Lecture 7

Database Management Systems

Winter 2004

CMPUT 391: Implementing Isolation

Dr. Osmar R. Zaïane

University of Alberta

Chapter 23 of Textbook

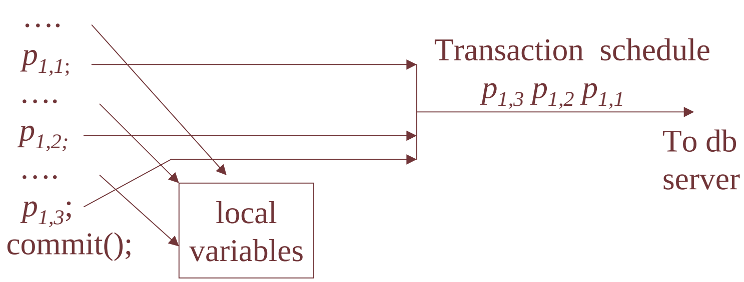
Based on slides by Lewis, Bernstein and Kifer.

Isolation

- Serial execution:
  - Since each transaction is consistent and isolated from all others, schedule is guaranteed to be correct for all applications
  - Inadequate performance
    - Since system has multiple asynchronous resources and transaction uses only one at a time
- Concurrent execution:
  - Improved performance (multiprogramming)
  - Some interleavings produce incorrect result
  - We are interested in concurrent schedules that are *equivalent* to serial schedules. These are referred to as *serializable* schedules.

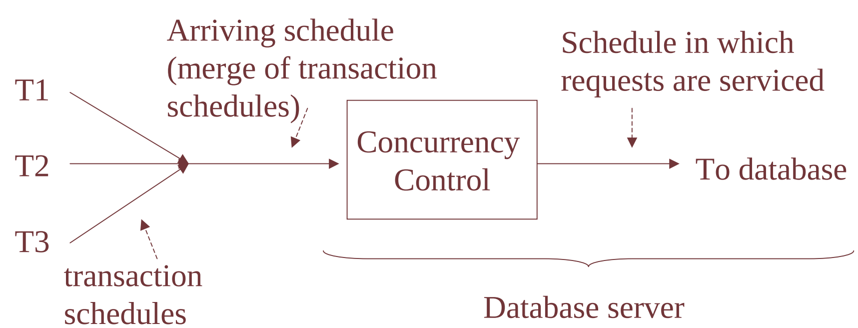
Transaction Schedule

T1: begin_transaction();



- Consistent - performs correctly when executed in isolation starting in a consistent database state
  - Preserves database consistency
  - Moves database to a new state that corresponds to new real-world state

Schedule



Schedule

- Representation 1:

T$_1$: p_1 p_2 p_3 p_4
T$_2$: p_1 p_2

time →

- Representation 2:

$p_{1,1}$ $p_{1,2}$ $p_{2,1}$ $p_{1,3}$ $p_{2,2}$ $p_{1,4}$
time →

Concurrency Control

- Transforms arriving schedule into a correct interleaved schedule to be submitted to the DBMS
  - Delays servicing a request (reordering) - causes a transaction to wait
  - Refuses to service a request - causes transaction to abort
- Actions taken by concurrency control have performance costs
  - Goal is to avoid delaying servicing a request

The Inconsistent Analysis Problem

- Occurs when a transaction reads several values from a database while a second transaction updates some of them.

T1	T2	A	B	C	sum
sum=0		$100	$50	$25	0
R(A)	R(A)	$100	$50	$25	0
sum=sum+A	A=A-10	$100	$50	$25	100
R(B)	W(A)	$90	$50	$25	100
sum=sum+B	R(C)	$90	$50	$25	150
	C=C+10	$90	$50	$25	150
	W(C)	$90	$50	$35	150
R(C)		$90	$50	$35	150
sum=sum+C		$90	$50	$35	185

Should be 175

Correct Schedules

- Interleaved schedules equivalent to serial schedules are the only ones guaranteed to be correct for *all* applications
- Equivalence based on *commutativity* of operations
- **Definition:** Database operations p_1 and p_2 commute if, for all initial database states, they return the same results and leave the database in the same final state when executed in either order.

Commutativity of Conventional Operations

- Read
  - *r(x, X)* - copy the value of database variable *x* to local variable *X*
- Write
  - *w(x, X)* - copy the value of local variable *X* to database variable *x*
- We use $r_1(x)$ and $w_1(x)$ to mean a read or write of *x* by transaction T_1

Commutativity of Read and Write Operations

- p_1 commutes with p_2 if
  - They operate on different data items
    - $w_1(x)$ commutes with $w_2(y)$ and $r_2(y)$
  - Both are reads
    - $r_1(x)$ commutes with $r_2(x)$
- Operations that do not commute *conflict*
  - $w_1(x)$ conflicts with $w_2(x)$
  - $w_1(x)$ conflicts with $r_2(x)$

	Read(x)	Write(x)
Read(x)	No	Yes
Write(x)	Yes	Yes

Equivalence of Schedules

- An interchange of adjacent operations of different transactions in a schedule creates an equivalent schedule if the operations commute

  $S_1 : S_{1,1}, p_{i,j}, p_{k,l}, S_{1,2}$ where $i \neq k$

  $S_2 : S_{1,1}, p_{k,l}, p_{i,j}, S_{1,2}$
- Equivalence is transitive: If S_1 is equivalent to S_2 (by a series of such interchanges), and S_2 is equivalent to S_3, then S_1 is equivalent to S_3

Example of Equivalence

conflict

S_1: $r_1(x)$ $r_2(x)$ $w_2(x)$ $r_1(y)$ $w_1(y)$

S_2: $r_1(x)$ $r_2(x)$ $r_1(y)$ $w_2(x)$ $w_1(y)$

S_3: $r_1(x)$ $r_1(y)$ $r_2(x)$ $w_2(x)$ $w_1(y)$

S_4: $r_1(x)$ $r_1(y)$ $r_2(x)$ $w_1(y)$ $w_2(x)$

S_5: $r_1(x)$ $r_1(y)$ $w_1(y)$ $r_2(x)$ $w_2(x)$

conflicting operations ordered in same way

S_1 is equivalent to S_5

S_5 is the serial schedule T_1, T_2

S_1 is *serializable*

S_1 is *not* equivalent to the serial schedule T_2, T_1

Example of Equivalence

T_1: *begin transaction*
read (x, X);
X = X+4;
write (x, X);
commit;

T_2: *begin transaction*
read (x,Y);
write (y,Y);
commit;

Interchange commuting operations

$x=1, y=3$ — $r_1(x)\ r_2(x)\ w_2(y)\ w_1(x)$ — $x=5, y=1$

$r_2(x)\ w_2(y)\ r_1(x)\ w_1(x)$ — $x=5, y=1$

T_2 T_1

Interchange conflicting operations

$x=1, y=3$ — $r_1(x)\ r_2(x)\ w_2(y)\ w_1(x)$ — $x=5, y=1$

$r_1(x)\ w_1(x)\ r_2(x)\ w_2(y)$ — $x=5, y=5$

T_1 T_2

Serializable Schedules

- S is serializable if it is equivalent to a serial schedule
- Transactions are totally isolated in a serializable schedule
- A schedule is correct for *any* application if it is a serializable schedule of consistent transactions
- The schedule :
  $r_1(x)\ r_2(y)\ w_2(x)\ w_1(y)$
  is *not* serializable

Isolation Levels

- Serializability provides a *conservative* definition of correctness
  - For a particular application there might be many acceptable *non*-serializable schedules
  - Requiring serializability might degrade performance
- DBMSs offer a variety of isolation levels:
  - SERIALIZABLE is the most stringent
  - Lower levels of isolation give better performance
    - *Might* allow incorrect schedules
    - *Might* be adequate for some applications

Serializable

- **Theorem -** Schedule S_1 can be derived from S_2 by a sequence of commutative interchanges if and only if conflicting operations in S_1 and S_2 are ordered in the same way

  If: A sequence of commutative interchanges can be determined that takes S_1 to S_2 since conflicting operations do not have to be reordered

  Only if: Commutative interchanges do not reorder conflicting operations

Conflict Equivalence

- **Definition**- Two schedules, S_1 and S_2, of the same set of operations are *conflict equivalent* if conflicting operations are ordered in the same way in both
  – Or (using theorem) if one can be obtained from the other by a series of commutative interchanges

Conflict Equivalence

- **Result**- A schedule is serializable if it is conflict equivalent to a serial schedule

$$r_1(x)\ w_2(x)\ w_1(y)\ r_2(y) \rightarrow r_1(x)\ w_1(y)\ w_2(x)\ r_2(y)$$

conflict ($r_1(x)$, $w_2(x)$) *conflict* ($w_1(y)$, $r_2(y)$)

- If in S transactions T_1 and T_2 have several pairs of conflicting operations ($p_{1,1}$ conflicts with $p_{2,1}$ and $p_{1,2}$ conflicts with $p_{2,2}$) then $p_{1,1}$ must precede $p_{2,1}$ and $p_{1,2}$ must precede $p_{2,2}$ (or vice versa) in order for S to be serializable.

Conflict Equivalence and Serializability

- Serializability is a conservative notion of correctness and conflict equivalence provides a conservative technique for determining serializability
- However, a concurrency control that guarantees conflict equivalence to serial schedules ensures correctness and is easily implemented

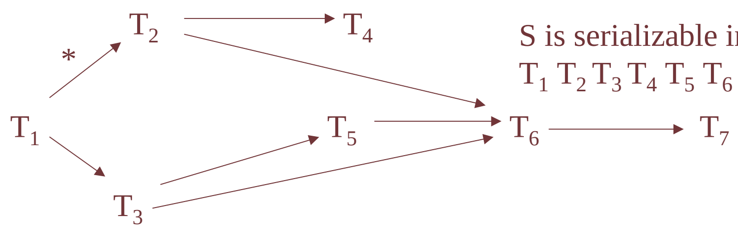
Serialization Graph of a Schedule, S

- Nodes represent transactions
- There is a directed edge from node T_i to node T_j if T_i has an operation $p_{i,k}$ that conflicts with an operation $p_{j,r}$ of T_j and $p_{i,k}$ precedes $p_{j,r}$ in S
- **Theorem** - A schedule is conflict serializable if and only if its serialization graph has no cycles

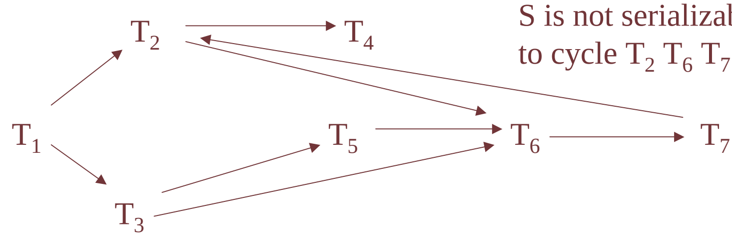
Example

Conflict (*)

S: … $p_{1,i}$, …, $p_{2,j}$, ...



S is serializable in order $T_1\ T_2\ T_3\ T_4\ T_5\ T_6\ T_7$

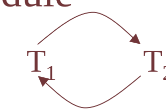


S is not serializable due to cycle $T_2\ T_6\ T_7\ T_2$

Intuition: Serializability and Nonserializability

- Consider the nonserializable schedule

$$r_1(x)\ w_2(x)\ r_2(y)\ w_1(y)$$



- Two ways to think about it:
  - Because of the read and write conflicts, the operations of T_1 and T_2 cannot be interchanged to make an equivalent serial schedule
  - Because T_1 read *x* before T_2 wrote it, T_1 must precede T_2 in any ordering, and because T_1 wrote *y* after T_2 read it, T_1 must follow T_2 in any ordering --- clearly an impossibility

Recoverability: Schedules with Aborted Transactions

T_1 : *r (x) w(y) commit*
T_2: *w(x)* *abort*

- T_2 has aborted but has had an indirect effect on the database – schedule is *un*recoverable
- **Problem**: T_1 read uncommitted data - *dirty read*
- **Solution:** A concurrency control is *recoverable* if it does not allow T_1 to commit until all other transactions that wrote values T_1 read have committed

T_1 : *r (x) w(y) req_commit* *abort*
T_2: *w(x)* *abort*

Cascaded Abort

- Recoverable schedules solve abort problem but allow *cascaded abort*: abort of one transaction forces abort of another

T_1: *r (y) w(z)* *abort*
T_2: *r (x) w(y)* *abort*
T_3: *w(x)* *abort*

- Better solution: prohibit dirty reads

Dirty Write

- *Dirty write*: A transaction writes a data item written by an active transaction
- Dirty write complicates rollback:

no rollback necessary ↓

T_1: *w(x)* *abort*
T_2 : *w(x)* *abort*

↑ what value of x should be restored?

Strict Schedules

- *Strict schedule*: Dirty writes and dirty reads are prohibited
- Strict and serializable are two different properties
  - Strict, non-serializable schedule:
    $r_1(x)\ w_2(x)\ r_2(y)\ w_1(y)\ c_1\ c_2$
  - Serializable, non-strict schedule:
    $w_2(x)\ r_1(x)\ w_2(y)\ r_1(y)\ c_1\ c_2$

Concurrency Control

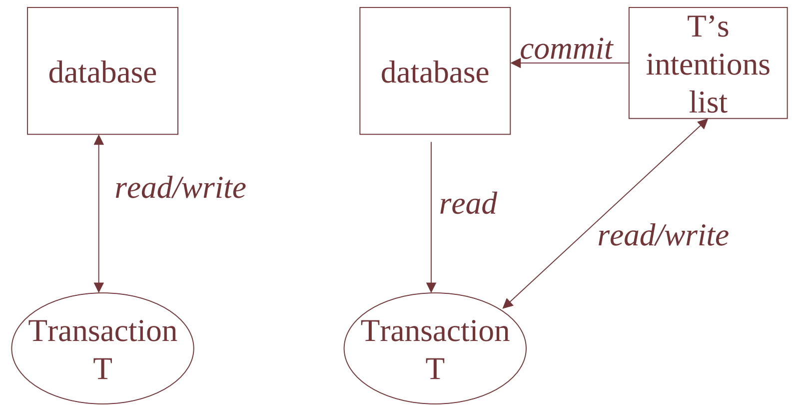
Arriving schedule (from transactions) → Concurrency Control → Serializable schedule (to processing engine)

- Concurrency control cannot see entire schedule:
  - It sees one request at a time and must decide whether to allow it to be serviced
- Strategy: Do not service a request if:
  - It violates strictness or serializability, or
  - There is a possibility that a subsequent arrival might cause a violation of serializability

Models of Concurrency Controls

- **Immediate Update**
  - A write updates a database item
  - A read copies value from a database item
  - Commit makes updates durable
  - Abort undoes updates
- **Deferred Update** – ***(we will likely not discuss this)***
  - A write stores new value in the transaction's intentions list (does not update database)
  - A read copies value from database or transaction's intentions list
  - Commit uses intentions list to durably update database
  - Abort discards intentions list

Immediate vs. Deferred Update



Immediate Update   Deferred Update

Models of Concurrency Controls

- **Pessimistic** –
  - A transaction requests permission for each database (read/write) operation
  - Concurrency control can:
    - *Grant* the operation (submit it for execution)
    - *Delay* it until a subsequent event occurs (commit or abort of another transaction), or
    - *Abort* the transaction
  - Decisions are made *conservatively* so that a commit request can *always* be granted
    - Takes precautions even if conflicts do not occur

Models of Concurrency Controls

- **Optimistic** -
  - Request for database operations (read/write) are *always* granted
  - Request to commit *might be denied*
    - Transaction is aborted if it performed a non-serializable operation
    - Assumes that conflicts are not likely
  - The earlier it can aborted the better

Deadlock

- **Problem**: Controls that cause transactions to wait can cause deadlocks

  $w_1(x)\ w_2(y)\ \ request\_r_1(y)\ \ request\_r_2(x)$
- **Solution**: Abort a transaction in the cycle
  - Use wait-for graph to detect cycle when a request is delayed or
  - Assume a deadlock when a transaction waits longer than some time-out period

Deadlock Prevention

- Assign priorities based on timestamps (i.e. The oldest transaction has higher priority).
- Assume T_i wants a lock that T_j holds. Two policies are possible:
  - Wait-Die: If T_i has higher priority, T_i allowed to wait for T_j; otherwise (T_i younger) T_i aborts
  - Wound-wait: If T_i has higher priority, T_j aborts; otherwise (T_i younger) T_i waits
- If a transaction re-starts, make sure it has its original timestamp

Deadlock and Timeouts

- A simple approach to deadlock prevention (and pseudo detection) is based on lock timeouts
- After requesting a lock on a locked data object, a transaction waits, but if the lock is not granted within a period (timeout), a deadlock is assumed and the waiting transaction is aborted and re-started.
- Very simple practical solution adopted by many DBMSs.

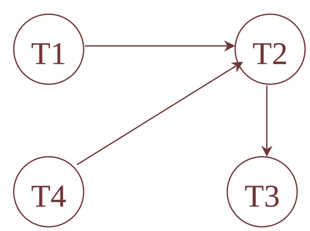
Deadlock Detection

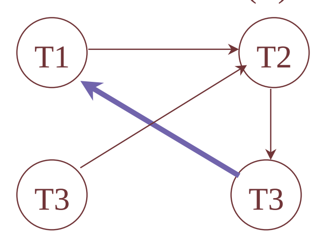
- Create a waits-for graph:
  - Nodes are transactions
  - There is an edge from T_i to T_j if T_i is waiting for T_j to release a lock
- Deadlock exists if there is a cycle in the graph.
- Periodically check for cycles in the waits-for graph.

Deadlock Detection (Continued)

Example:

T1:	S(A), R(A),		S(B)				
T2:		X(B),W(B)			X(C)		
T3:				S(C), R(C)			X(A)
T4:						X(B)	





Locking Implementation of an Immediate-Update Pessimistic Control

- A transaction can read a database item if it holds a read (shared) lock on the item
- It can read *or* update the item if it holds a write (exclusive) lock
- If the transaction does not already hold the required lock, a lock request is automatically made as part of the access

Locking

- Request for read lock granted if no transaction currently holds write lock on item
  - Cannot read an item written by an active transaction
- Request for write lock granted if no transaction holds any lock on item
  - Cannot write an item read/written by an active transaction

	Granted mode	
Requested mode	*read*	*write*
read		x
write	x	x

Locking

- All locks held by a transaction are released when the transaction completes (commits or aborts)

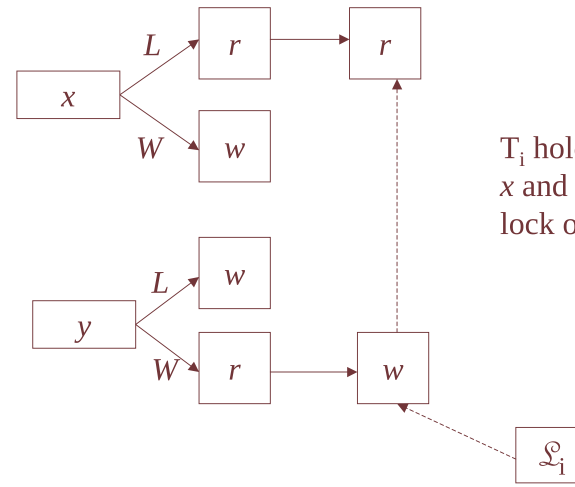
Locking

- **Result**: A lock is not granted if the requested access conflicts with a prior access of an active transaction; instead the transaction waits. This enforces the rule:
  - Do not grant a request that imposes an ordering among active transactions (delay the requesting transaction)
- Resulting schedules are serializable and strict

Locking Implementation

- Associate a *lock set, L(x)*, and a *wait set, W(x)*, with each active database item, *x*
  - *L(x)* contains an entry for each granted lock
  - *W(x)* contains an entry for each pending request
  - When an entry is removed from *L(x)* (due to transaction termination), promote (non-conflicting) entries from *W(x)* using some scheduling policy (*e.g.*, FCFS)
- Associate a lock list, $\mathcal{L}_i$, with each transaction, T_i.
  - $\mathcal{L}_i$ links T_i's elements in all lock and wait sets
  - Used to release locks on termination

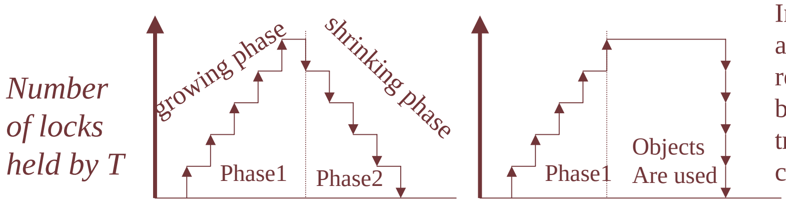
Locking Implementation



T_i holds an *r* lock on *x* and waits for a *w* lock on *y*

Two-Phase Locking

- Transaction does not release a lock until it has all the locks it will ever require.
- Transaction, T, has a locking phase followed by an unlocking phase



2-phase locking (2PL) Strict 2-phase locking (strict 2PL)

In strict-2PL all locks are released at one before the transaction commits

- Guarantees serializability when locking is done manually

Two-Phase Locking

- **Theorem**: A concurrency control that uses two phase locking produces only serializable schedules.
  - *Proof:* Consider two transactions T_1 and T_2 in schedule S produced by a two-phase locking control and assume T_1's first unlock precedes T_2's first unlock.
    - If they do not access common data items, then all operations commute and S is serializable.
    - Suppose they do. For each common item *x*, all of T_1's accesses to *x* precede all of T_2's. If this were not the case, T_2's first unlock must precede a lock request of T_1. Since both transactions are two-phase, this implies that T_2's first unlock precedes T_1's first unlock, contradicting the assumption.
    - Thus S is serializable.

Two-Phase Locking

- A schedule produced by a two-phase locking control is:
  - Equivalent to a serial schedule in which transactions are ordered by the time of their first unlock operation
  - Not necessarily recoverable (dirty reads and writes are possible)

T1: *l(x) r(x) l(y) w(y) u(y)* *abort*
T2: *l(y) r(y) l(z) w(z) u(z) u(y) commit*

Two-Phase Locking

- A two-phase locking control that holds write locks until commit produces strict serializable schedules
- A strict two-phase locking control holds all locks until commit and produces strict serializable schedules
  - This is automatic locking
  - Equivalent to a serial schedule in which transactions are ordered by their commit time
- "Strict" is used in two different ways: a control that releases read locks early guarantees *strictness*, but is not *strict* two-phase locking control

Lock Granularity

- Data item: variable, record, row, table, file
- When an item is accessed, the DBMS locks an entity that contains the item. The size of that entity determines the *granularity* of the lock
  - Coarse granularity (large entities locked)
    - **Advantage**: If transactions tend to access multiple items in the same entity, fewer lock requests need to be processed and less lock storage space required
    - **Disadvantage:** Concurrency is reduced since some items are unnecessarily locked
  - Fine granularity (small entities locked)
    - Advantages and disadvantages are reversed

Lock Granularity

- Table locking (*coarse*)
  - Lock entire table when a row is accessed.
- Row (tuple) locking (*fine*)
  - Lock only the row that is accessed.
- Page locking (compromise)
  - When a row is accessed, lock the containing page

Timestamp-Ordered Concurrency Control

- Each transaction given a (unique) timestamp (current clock value) when initiated
- Uses the immediate update model
- Guarantees equivalent serial order based on timestamps (initiation order)
  - Control is *static* (as opposed to *dynamic*, in which the equivalent serial order is determined as the schedule progresses)

Timestamp-Ordered Concurrency Control

- Associated with each database item, *x*, are two timestamps:
  - *wt(x)*, the largest timestamp of any transaction that has written *x*,
  - *rt(x)*, the largest timestamp of any transaction that has read *x*,
  - and an indication of whether or not the last write to that item is from a committed transaction

Timestamp-Ordered Concurrency Control

- If T requests to read *x*:
  - **R1**: if $TS(T) < wt(x)$, then T is too old; abort T
  - **R2**: if $TS(T) > wt(x)$, then
    - if the value of *x* is committed, grant T's read and if $TS(T) > rt(x)$ assign $TS(T)$ to $rt(x)$
    - if the value of *x* is not committed, T waits (to avoid a dirty read)

Timestamp-Ordered Concurrency Control

- If T requests to write *x* :
  - **W1**: If $TS(T) < rt(x)$, then T is too old; abort T
  - **W2**: If $rt(x) < TS(T) < wt(x)$, then no transaction that read *x* should have read the value T is attempting to write and no transaction will read that value (R1)
    - If *x* is committed, grant the request but do not do the write
    - If *x* is not committed, T waits to see if newer value will commit. If it does, discard T's write, else perform it
  - **W3**: If $wt(x), rt(x) < TS(T)$, then if *x* is committed, grant the request and assign $TS(T)$ to $wt(x)$, else T waits

Example

- Assume $TS(T_1) < TS(T_2)$, at t_0 x and y are committed, and x's and y's read and write timestamps are less than $TS(T_1)$

T_1 :		*r(y)*				*w(x)*	*commit*
T_2:			*w(y)*	*w(x)*	*commit*		
	t_0	t_1	t_2	t_3		t_4	

t_1: (R2) $TS(T_1) > wt(y)$; assign $TS(T_1)$ to $rt(y)$

t_2: (W3) $TS(T2) > rt(y), wt(y)$; assign $TS(T_2)$ to $wt(y)$

t_3: (W3) $TS(T2) > rt(x), wt(x)$; assign $TS(T_2)$ to $wt(x)$

t_4: (W2) $rt(x) < TS(T1) < wt(x)$; grant request, but don't do the write

Timestamp-Ordered Concurrency Control

- Control accepts schedules that are *not conflict equivalent* to any serial schedule and would not be accepted by a two-phase locking control
  - Previous example equivalent to T_1, T_2
- But additional space required in database for storing timestamps and time for managing timestamps
  - Reading a data item now implies writing back a new value of its timestamp

Optimistic Concurrency Control

- No locking (and hence no waiting) means deadlocks are not possible
- Rollback is a problem if optimistic assumption is not valid: work of entire transaction is lost
  - With two-phase locking, rollback occurs only with deadlock
  - With timestamp-ordered control, rollback is detected before transaction completes